# A Local Computation Approximation Scheme to Maximum Matching 

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#### Abstract

We present a polylogarithmic local computation matching algorithm which guarantees a (1- $\epsilon$ )-approximation to the maximum matching in graphs of bounded degree.


## 1 Introduction

Finding matchings - sets of vertex disjoint edges in a graph - has been an important topic of research for computer scientists for over 50 years. Of particular importance is finding maximum matchings - matchings of maximal cardinality. Algorithms that find a maximum matching have many applications in computer science; in fact, their usefulness extends far beyond the boundaries of computer science - to disciplines such as economics, biology and chemistry.

The first works on matching were based on unweighted bipartite graphs (representing problems such as matching men and women). Hall's marriage theorem [6] gives a necessary and sufficient condition for the existence of a perfect matching ${ }^{1}$. The efficient algorithms for the weighted bipartite matching problem date back to the Hungarian method [12|18]. In this work we focus on maximum matchings in general unweighed graphs. Berge [3] proved that a matching is a maximum matching if and only if the graph has no augmenting paths with respect to the matching. Edmonds used augmenting paths to find a maximum matching in his seminal work [5], in which he showed that a maximum matching can be found in polynomial time. Much work on matching been done since (e.g., [7/9]16[17]). Our work uses ideas from Hopcroft and Karp's algorithm for finding maximal matching in bipartite graphs [9], which runs in time $O\left(n^{2.5}\right)$.

Local computation algorithms (LCAs) [20] consider the scenario in which we must respond to queries (regarding a feasible solution) quickly and efficiently, yet we never need the entire solution at once. The replies to the queries need to be consistent; namely, the responses to all possibly queries combine to a single feasible solution. For example, an LCA for matching in a graph $G$, receives an edge-query for an edge $e \in G$ and replies "yes" if and only if $e$ is part of the matching. The replies to all the possible edge queries define a matching in the graph.

In this work we present a local computation approximation scheme to maximum matching. Specifically, we present an LCA such that for any $\epsilon>0$, the edge-query replies comprise a matching that is a $(1-\epsilon)$-approximation to the maximum matching. Our LCA requires $O\left(\log ^{3} n\right)$ space, and with probability at least $1-1 / n^{2}$, for any edge-query, it runs in time $O\left(\log ^{4} n\right)$. To the best of our knowledge, this is the first local computation approximation algorithm for a problem which provably does not have an LCA.

Related work. In the distributed setting, Itai and Israeli [10] showed a randomized algorithm which computes a maximal matching (which is a $1 / 2$-approximation to the maximum matching) and runs in $O(\log n)$ time with high probability. This result has been improved several times since (e.g., [4]8]); of particular relevance is the approximation scheme of Lotker et al. [13], which, for every $\epsilon>0$, computes a $(1-\epsilon)$-approximation to the maximum matching in $O(\log n)$ time. Kuhn et al., [11] proved that any distributed algorithm, randomized or deterministic, requires (in expectation) $\Omega(\sqrt{\log n / \log \log n})$ time to compute a $\Theta(1)$-approximation to the maximum matching, even if the message size is unbounded.

[^0]Rubinfeld et al., [20] showed how to transform distributed algorithms to LCAs, and gave LCAs for several problems, including maximal independent set and hypergraph 2-coloring. Unfortunately, their method bounds the running time of the transformed algorithm exponentially in the running time of the distributed algorithm. Therefore, distributed algorithms for approximate maximum matching cannot be (trivially) transformed to LCAs using their technique.

Query trees model the dependency of queries on the replies to other queries, and were introduced in the local setting by Nguyen and Onak [19]. If a random permutation of the vertices is generated, and a sequential algorithm is simulated on this order, the reply to a query on vertex $v$ depends only on the replies to queries on the neighbors of $v$ which come before it in the permutation. Alon et al., [2] showed that if the running time of an algorithm is $O(f(n))$, where $f$ is polylogarithmic in $n$, a $1 / n^{2}$ - almost $f(n)$-independent ordering on the vertices can be generated in time $O\left(f(n) \log ^{2} n\right)$, thus guaranteeing the polylogarithmic space bound of any such algorithm. Mansour et al., [14] showed that the size of the query tree can be bounded, with high probability, by $O(\log n)$, for graphs of bounded degree. They also showed that it is possible to transform many on-line algorithms to LCAs. One of their examples is an LCA for maximal matching, which immediately gives a $1 / 2$-approximation to the maximum matching. In a recent work, [15], LCAs were presented for mechanism design problems. One of their impossibility results shows that any LCA for maximum matching requires $\Omega(n)$ time.

## 2 Notation and Preliminaries

### 2.1 Graph Theory

For an undirected graph $G=(V, E)$, a matching is a subset of edges $M \subseteq E$ such that no two edges $e_{1}, e_{2} \in M$ share a vertex. We denote by $M^{*}$ a matching of maximum cardinality. An augmenting path with respect to a matching $M$ is a simple path whose endpoints are free (i.e., not part of any edge in the matching $M$ ), and whose edges alternate between $E \backslash M$ and $M$. A set of augmenting paths $P$ is independent if no two paths $p_{1}, p_{2} \in P$ share a vertex.

For sets $A$ and $B$, we denote $A \oplus B \stackrel{\text { def }}{=}(A \cup B) \backslash(A \cap B)$. An important observation regarding augmenting paths and matchings is the following.

Observation 1 If $M$ is a matching and $P$ is an independent set of augmenting paths, then $M \oplus P$ is a matching of size $|M|+|P|$.

A vertex $u \in V$ is a neighbor of vertex $v \in V$ if $(u, v) \in E$. Let $N(v)$ denote the set of neighbors of $v$, i.e., $N(v)=\{u:(v, u) \in E\}$. We assume that we have direct access both to $N(v)$ and to individual edges.

An independent set (IS) is a subset of vertices $W \subseteq V$ with the property that for any $u, v \in W$ we have $(u, v) \notin E$, namely, no two vertices $u, v \in W$ are neighbors in $G$. The IS is maximal (denoted by MIS) if no other vertices can be added to it without violating the independence property.

### 2.2 Local Computation Algorithms

We use the following model of local computation algorithms (LCAs) $20{ }^{2} \mathrm{~A}(t(n), s(n), \delta(n))$ - local computation algorithm $\mathcal{L} \mathcal{A}$ for a computational problem is a (randomized) algorithm which receives an input of size $n$, and a query $x$. Algorithm $\mathcal{L} \mathcal{A}$ uses at most $s(n)$ memory, and with probability at least $1-\delta(n)$, it replies to any query $x$ in time $t(n)$. The algorithm must be consistent, that is, the replies to all of the possible queries combine to a single feasible solution to the problem.

[^1]
### 2.3 Query Trees

Let $G=(V, E)$ be a graph of bounded degree $d$. A real number $r(v) \in[0,1]$ is assigned independently and uniformly at random to every vertex $v$ in the graph. We refer to this random number as the rank of $v$. Each vertex in the graph $G$ holds an input $x(v) \in R$, where the range $R$ is some finite set. A randomized Boolean function $F$ is defined inductively on the vertices in the graph such that $F(v)$ is a function of the input $x(v)$ at $v$ as well as the values of $F$ at the neighbors $w$ of $v$ for which $r(w)<r(v)$.

We would like to upper bound the number of queries that are needed to be made vertices in the graph in order to compute $F\left(v_{0}\right)$ for any vertex $v_{0} \in G$. We turn to the simpler task of bounding the size of a certain $d$-regular tree, which is an upper bound on the number of queries. Consider an infinite $d$-regular tree $\mathcal{T}$ rooted at $v_{0}$. Each node $w$ in $\mathcal{T}$ is assigned independently and uniformly at random a distinct real number $r(w) \in[0,1]$. For every node $w \in \mathcal{T}$ other than $v_{0}$, let parent $(w)$ denote the parent node of $w$. We grow a (possibly infinite) subtree $T$ of $\mathcal{T}$ rooted at $v$ as follows: a node $w$ is in the subtree $T$ if and only if parent $(w)$ is in $T$ and $r(w)<r($ parent $(w))$. We keep growing $T$ in this manner such that a node $w^{\prime} \in T$ is a leaf node in $T$ if the ranks of its $d$ children are all larger than $r\left(w^{\prime}\right)$. We call the random tree $T$ constructed in this way a query tree and we denote by $|T|$ the random variable that corresponds to the size of $T$. Note that $|T|$ is an upper bound on the number of queries.

If the reply to a query $q$ depends (only) on the replies to a set of queries, $Q$, we call $Q$ the set of relevant queries with respect to $q$.

### 2.4 Random Orders

Let $[n]$ denote the set $\{1, \ldots, n\}$.
A distribution $D:\{0,1\}^{n} \rightarrow \mathbb{R}^{\geq 0}$ is $k$-wise independent if, when $D$ is restricted to any index subset $S \subset[n]$ of size at most $k$, the induced distribution over $S$ is the uniform distribution.

A random ordering $D_{\mathbf{r}}$ induces a probability distribution over permutations of $[n]$. It is said to $\epsilon$-almost $k$-wise independent if for any subset $S \subset[n]$ of size at most $k$, the variation distance between the distribution induced by $D_{\mathbf{r}}$ on $S$ and a uniform permutation over $S$ is at most $\epsilon$. We use the following Theorem from [2].

Theorem 2 ([|2]). Let $n \geq 2$ be an integer and let $2 \leq k \leq n$. Then there is a construction of $\frac{1}{n^{2}}$-almost $k$-wise independent random ordering over $[n]$ whose seed length is $O\left(k \log ^{2} n\right)$.

We provide a short, intuitive explanation of the construction. We can construct $n k$-wise independent random variables $Z=\left(z_{1}, \ldots, z_{n}\right)$, using a seed of length $k \log n$ (see [1]). We generate $4 \log n$ independent copies of $k$-wise independent random variables, $Z_{1}, \ldots Z_{4} \log n$. For $i \in[n]$, taking the $i$-th bit of each $Z_{j}, 1 \leq j \leq 4 \log n$ makes for a random variable $r(i) \in\{0,1\}^{4 \log n}$, which can be expressed as an integer in $\left\{0,1, \ldots, n^{4}-1\right\}$. The order is induced by $r$ ( $u$ comes before $v$ in the order if $r(u)<r(v)$ ). The probability that there exists $u, v \in[n]$ such that $r(u)=r(v)$ is at most $1 / n^{2}$, hence the ordering is $1 / n^{2}$-almost $k$-wise independent.

## 3 Approximate Maximum Matching

We present a local computation approximation scheme for maximum matching: We show an LCA that, for any $\epsilon>0$, computes a maximal matching which is a $(1-\epsilon)$-approximation to the maximum matching.

Our main result is the following theorem:
Theorem 3. Let $G=(V, E)$ be a graph of bounded degree d. Then there exists an $\left(O\left(\log ^{4} n\right), O\left(\log ^{3} n\right), 1 / n\right)$ LCA that, for every $\epsilon>0$, computes a maximal matching which is a $(1-\epsilon)$-approximation to the maximum matching.

Our algorithm is, in essence, an implementation of the abstract algorithm of Lotker et al., [13]. Their algorithm, relies on several interesting results due to Hopcroft and Karp [9]. First, we briefly recount some of these results, as they are essential for the understanding of our algorithm.

### 3.1 Distributed Maximal Matching

While the main result of Hopcroft and Karp [9] is an improved matching algorithm for bipartite graphs, they show the following useful lemmas for general graphs. The first lemma shows that if the current matching has augmenting paths of length at least $\ell$, then using a maximal set of augmenting paths of length $\ell$ will result in a matching for which the shortest augmenting path is strictly longer than $\ell$. This gives a natural progression for the algorithm.

Lemma 4. [9] Let $G=(V, E)$ be an undirected graph, and let $M$ be some matching in $G$. If the shortest augmenting path with respect to $M$ has length $\ell$ and $\Phi$ is a maximal set of independent augmenting paths of length $\ell$, the shortest augmenting path with respect to $M \oplus \Phi$ has length strictly greater than $\ell$.

The second lemma shows that if there are no short augmenting paths then the current matching is approximately optimal.

Lemma 5. [9] Let $G=(V, E)$ be an undirected graph. Let $M$ be some matching in $G$, and let $M^{*}$ be a maximum matching in $G$. If the shortest augmenting path with respect to $M$ has length $2 k-1>1$ then $|M| \geq(1-1 / k)\left|M^{*}\right|$.

Lotker et al., [13] gave the following abstract approximation scheme for maximal matching in the distributed setting ${ }^{3}$ Start with an empty matching. In stage $\ell=1,3, \ldots, 2 k-1$, add a maximal independent collection of augmenting paths of length $\ell$. For $k=\lceil 1 / \epsilon\rceil$, by Lemma[5] we have that the matching $M_{\ell}$ is a $(1-\epsilon)$-approximation to the maximum matching.

In order to find such a collection of augmenting paths of length $\ell$, we need to define a conflict graph:
Definition 6. [13] Let $G=(V, E)$ be an undirected graph, let $M \subseteq E$ be a matching, and let $\ell>0$ be an integer. The $\ell$-conflict graph with respect to $M$ in $G$, denoted $C_{M}(\ell)$, is defined as follows. The nodes of $C_{M}(\ell)$ are all augmenting paths of length $\ell$, with respect to $M$, and two nodes in $C_{M}(\ell)$ are connected by an edge if and only if their corresponding augmenting paths intersect at a vertex of $G \square$

We present the abstract distributed algorithm of [13], AbstractDistributedMM.

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Algorithm 1 - AbstractDistributedMM - Abstract distributed algorithm with input \(G=(V, E)\) and \(\epsilon>0\)
    \(M_{-1} \leftarrow \emptyset\)
    \(k \leftarrow\lceil 1 / \epsilon\rceil\)
    for \(\ell \leftarrow 1,3, \ldots, 2 k-1\), do
        Construct the conflict graph \(C_{M_{\ell-2}}(\ell)\)
        Let \(\mathcal{I}\) be an MIS of \(C_{M_{\ell-2}}(\ell)\)
        Let \(\Phi\left(M_{\ell-2}\right)\) be the union of augmenting paths corresponding to \(\mathcal{I}\)
        \(M_{\ell} \leftarrow M_{\ell-2} \oplus \Phi\left(M_{\ell-2}\right) \quad \triangleright M_{\ell}\) is matching at the end of phase \(\ell\)
    end for
    Output \(M_{\ell} \quad \triangleright M_{\ell}\) is a \(\left(1-\frac{1}{k+1}\right)\)-approximate maximum matching
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Note that for $M_{\ell}$, the minimal augmenting path is of length at least $\ell+2$. This follows since $\Phi\left(M_{\ell-2}\right)$ is a maximal independent set of augmenting paths of length $\ell$. When we add $\Phi\left(M_{\ell-2}\right)$ to $M_{\ell-2}$, to get $M_{\ell}$, by Lemma 4 all the remaining augmenting paths are of length at least $\ell+2$ (recall that augmenting paths have odd lengths).

Lines 4 - 7 do the task of computing $M_{\ell}$ as follows: the conflict graph $C_{M_{\ell-2}}(\ell)$ is constructed and an MIS, $\Phi\left(M_{\ell-2}\right)$, is found in it. $\Phi\left(M_{\ell-2}\right)$ is then used to augment $M_{\ell-2}$, to give $M_{\ell}$.

We would like to simulate this algorithm locally. Our main challenge is to simulate Lines 4 - 7 without explicitly constructing the entire conflict graph $C_{M_{\ell-2}}(\ell)$. To do this, we will simulate an on-line MIS algorithm.

[^2]
### 3.2 Local Simulation of the On-Line Greedy MIS Algorithm

In the on-line setting, the vertices arrive in some unknown order, and GreedyMIS operates as follows: Initialize the set $I=\emptyset$. When a vertex $v$ arrives, GreedyMIS checks whether any of $v$ 's neighbors, $N(v)$, is in $I$. If none of them are, $v$ is added to $I$. Otherwise, $v$ is not in $I$. (The pseudocode for GreedyMIS can be found in the full version of the paper.)

In order to simulate GreedyMIS locally, we first need to fix the order (of arrival) of the vertices, $\pi$. If we know that each query depends on at most $k$ previous queries, we do not need to explicitly generate the order $\pi$ on all the vertices (as this would take at least linear time). By Theorem 2 we can produce a $\frac{1}{n^{2}}$-almost- $k$-wise independent random ordering on the edges, using a seed, $s$, of length $O\left(k \log ^{2} n\right)$.

Technically, this is done as follows. Let $r$ be a function $r:(v, s) \rightarrow\left[c n^{4}\right]$, for some constant $c{ }^{5}$ The vertex order $\pi$ is determined as follows: vertex $v$ appears before vertex $u$ in the order $\pi$ if $r(v, s)<r(u, s)$. Let $G^{\prime}=\left(V^{\prime}, E^{\prime}\right)$ be the subgraph of $G$ induced by the vertices $V^{\prime} \subseteq V$; we denote by $\pi\left(G^{\prime}, s\right)$ the partial order of $\pi$ on $V^{\prime}$. Note that we only need to store $s$ in the memory: we can then compute, for any subset $V^{\prime}$, the induced order of their arrival.

When simulating GreedyMIS on the conflict graph $C_{M}(\ell)=\left(V_{C_{M}}, E_{C_{M}}\right)$, we only need a subset of the nodes, $V^{\prime} \subseteq V_{C_{M}}$. Therefore, there is no need to construct $C_{M}(\ell)$ entirely; only the relevant subgraph need be constructed. This is the main observation which allows us to bound the space and time required by our algorithm.

### 3.3 LCA for Maximal Matching

We present our algorithm for maximal matching - LocalMM, and analyze it. (The pseudocode for LocalMM can be found in the full version of the paper.) In contrast to the distributed algorithm, which runs iteratively, LocalMM is recursive in nature. In each iteration of AbstractDistributedMM, a maximal matching $M_{\ell}$, is computed, where $M_{\ell}$ has no augmenting path of length less than $\ell$. We call each such iteration a phase, and there are a total of $k$ phases: $1,3, \ldots 2 k-1$. To find out whether an edge $e \in E$ is in $M_{\ell}$, we recursively compute whether it is in $M_{\ell-2}$ and whether it is in $\Phi\left(M_{\ell-2}\right)$, a maximal set of augmenting paths of length $\ell$. We use the following simple observation to determine whether $e \in M_{\ell}$. The observation follows since $M_{\ell} \leftarrow M_{\ell-2} \oplus \Phi\left(M_{\ell-2}\right)$.

Observation $7 e \in M_{\ell}$ if and only if it is in either in $M_{\ell-2}$ or in $\Phi\left(M_{\ell-2}\right)$, but not in both.
Recall that LocalMM receives an edge $e \in E$ as a query, and outputs "yes/no". To determine whether $e \in M_{2 k-1}$, it therefore suffices to determine, for $\ell=1,3, \ldots, 2 k-3$, whether $e \in M_{\ell}$ and whether $e \in \Phi\left(M_{\ell}\right)$.

We will outline our algorithm by tracking a single query. (The initialization parameters will be explained at the end.) When queried on an edge $e$, LocalMM calls the procedure ISInMATCHING with $e$ and the number of phases $k$. For clarity, we sometimes omit some of the parameters from the descriptions of the procedures.

Procedure ISInMATCHING determines whether an edge $e$ in in the matching $M_{\ell}$. To determine whether $e \in M_{\ell}$, IsInMatching recursively checks whether $e \in M_{\ell-2}$, by calling $\operatorname{IsInMATCHING}(\ell-2)$, and whether $e$ is in some path in the MIS $\Phi\left(M_{\ell-2}\right)$ of $C_{M_{\ell-2}}(\ell)$. This is done by generating all paths $p$ of length $\ell$ that include $e$, and calling IsPathinMIS $(p)$ on each. IsPathInMIS $(p)$ checks whether $p$ is an augmenting path, and if so, whether it in the independent set of augmenting paths. By Observation 7 we can compute whether $e$ is in $M_{\ell}$ given the output of the calls.

Procedure ISPATHINMIS receives a path $p$ and returns whether the path is in the MIS of augmenting paths of length $\ell$. The procedure first computes all the relevant augmenting paths (relative to $p$ ) using RELEVANTPaths. Given the set of relevant paths (represented by nodes) and the intersection between them (represented by edges) we simulate GreedyMIS on this subgraph. The resulting independent set is a set of independent augmenting paths. We then just need to check if the path $p$ is in that set.

[^3]Procedure RelevantPaths receives a path $p$ and returns all the relevant augmenting paths relative to $p$. The procedure returns the subgraph of $C_{M_{\ell-2}}(\ell), C=\left(V_{C}, E_{C}\right)$, which includes $p$ and all the relevant nodes. These are exactly the nodes needed for the simulation of GreedyMIS, given the order induced by seed $s_{\ell}$. The set of augmenting paths $V_{C}$ is constructed iteratively, by adding an augmenting path $q$ if it intersects some path $q^{\prime} \in V_{C}$ and arrives before it (i.e., $r\left(q, s_{\ell}\right)<r\left(q^{\prime}, s_{\ell}\right)$ ). In order to determine whether to add path $q$ to $V_{C}$, we need first to test if $q$ is indeed a valid augmenting path, which is done using ISANAUGMENTINGPATH.

Procedure IsAnAugmenting Path tests if a given path $p$ is an augmenting path. It is based on the following observation.

Observation 8 For any graph $G=(V, E)$, let $M$ be a matching in $G$, and let $p=e_{1}, e_{2}, \ldots, e_{\ell}$ be a path in $G$. Path $p$ is an augmenting path with respect to $M$ if and only if all odd numbered edges are not in $M$, all even numbered edges are in $M$, and both the vertices at the ends of $p$ are free.

Given a path $p$ of length $\ell$, to determine whether $p \in C_{M_{\ell-2}}(\ell)$, IsAnAUGMENTingPath $(\ell)$ determines, for each edge in the path, whether it is in $M_{\ell-2}$, by calling IsInMatching $(\ell-2)$. It also checks whether the end vertices are free, by calling Procedure $\operatorname{ISFreE}(\ell)$, which checks, for each vertex, if any of its adjacent edges are in $M_{\ell-2}$. From Observation 8, IsAnAUGMENTINGPath $(\ell)$ correctly determines whether $p$ is an augmenting with respect to $M_{\ell-2}$.

We end by describing the initialization procedure Initialize, which is run only once, during the first query. The procedure sets the number of phases to $\lceil 1 / \epsilon\rceil$. It is important to set a different seed $s_{\ell}$ for each phase $\ell$, since the conflict graphs are unrelated (and even the size of the description of each node, a path of length $\ell$, is different). The lengths of the $k$ seeds, $s_{1}, s_{3}, \ldots, s_{2 k-1}$, determine our memory requirement.

### 3.4 Bounding the Complexity

In this section we prove Theorem 3 We start with the following observation:
Observation 9 In any graph $G=(V, E)$ with bounded degree d, each edge $e \in E$ can be part of at most $\ell(d-1)^{\ell-1}$ paths of length $\ell$. Furthermore, given $e$, it takes at most $O\left(\ell(d-1)^{\ell-1}\right)$ time to find all such paths.

Proof. Consider a path $p=\left(e_{1}, e_{2}, \ldots, e_{\ell}\right)$ of length $\ell$. If $p$ includes the edge $e$, then $e$ can be in one of the $\ell$ positions. Given that $e_{i}=e$, there are at most $d-1$ possibilities for $e_{i+1}$ and for $e_{i-1}$, which implies at most $(d-1)^{\ell-1}$ possibilities to complete the path to be of length $\ell$.

Observation 9 yields the following corollary.
Corollary 10. The $\ell$-conflict graph with respect to any matching $M$ in $G=(V, E), C_{M}(\ell)$, consists of at most $\ell(d-1)^{\ell-1}|E|=O(|V|)$ nodes, and has maximal degree at most $d(\ell+1) \ell(d-1)^{\ell-1}$.

Proof. (For the degree bound.) Each path has length $\ell$, and therefore has $\ell+1$ vertices. Each vertex has degree at most $d$, which implies $d(\ell+1)$ edges. Each edge is in at most $\ell(d-1)^{\ell-1}$ paths.

Our main task will be to compute a bound on the number of recursive calls. First, let us summarize a recursive call. The only procedure whose runtime depends on the order induced by $s_{\ell}$ is RelevantPaths, which depends on the number of vertices $V_{C}$ (which is a random variable depending of the seed $s_{\ell}$ ). To simplify the notation we define the random variable $X_{\ell}=d(\ell+1) \ell(d-1)^{\ell-1}\left|V_{C}\right|$. Technically, GreedyMIS also depends on $V_{C}$, but its running time is dominated by the running time of RELEVANTPATHS.

| Calling procedure | Called Procedures |
| :---: | :---: |
| IsInMATCHING $(\ell)$ | $1 \times \operatorname{IsInMATCHING}(\ell-2)$ and $\ell(d-1)^{\ell-1} \times \operatorname{IsPATHINMIS}(\ell)$ |
| ISPATHINMIS $(\ell)$ | $1 \times \operatorname{RELEVANTPATHS}(\ell)$ and $1 \times \operatorname{GreedyMIS}$ |
| RELEVANTPATHS $(\ell)$ | $X_{\ell} \times \operatorname{ISANAUGMENTINGPATH}(\ell)$ |
| ISANAUGMENTINGPATH $(\ell)$ | $\ell \times \operatorname{IsInMATCHING}(\ell-2)$ and $2 \times \operatorname{IsFREE}(\ell)$ |
| ISFREE $(\ell)$ | $(d-1) \times \operatorname{ISInMATCHING}(\ell-2)$ |

From the table, it is easy to deduce the following proposition.

Proposition 11. IsAnAUGMENTINGPATH $(\ell)$ generates at most $\ell+2(d-1)$ calls to $\operatorname{IsInMATCHING}(\ell-2)$, and therefore at most $(\ell+2 d-2) \cdot \ell(d-1)^{\ell-1}$ calls to $\operatorname{ISPATHINMIS}(\ell-2)$.

We would like to bound $X_{\ell}$, the number of calls to IsAnAUGMENTINGPATH $(\ell)$ during a single execution of IsPathinMIS $(G, p, \ell, S)$. We require the following theorem, the proof of which appears in Section 4

Theorem 12. For any infinite query tree $T$ with bounded degree $d$, there exists a constant $c$, which depends only on $d$, such that for any large enough $N>0$,

$$
\operatorname{Pr}[|T|>N] \leq e^{-c N}
$$

As a query tree $T$ of bounded degree $D=d(\ell+1) \ell(d-1)^{\ell-1}$ is an upper bound to $X_{\ell}$ (by Corollary 10, $D$ is an upper bound on the degree of $C_{M_{\ell-2}}(\ell)$ ), we have the following corollary to Theorem 12 ,

Corollary 13. There exists an absolute constant $c$, which depends only on $d$, such that for any large enough $N>0$,

$$
\operatorname{Pr}\left[X_{\ell}>N\right] \leq e^{-c N}
$$

Denote by $f_{\ell}$ the number of calls to ISAnAUGMENTINGPATH $(\ell)$ during one execution of LocalMM. Let $f=$ $\sum_{\ell=1}^{2 k-1} f_{\ell}{ }^{6}$ The base cases of the recursive calls LocalMM makes are ISANAUGMENTINGPATH (1) (which always returns TRUE). As the execution of each procedure of LocalMM results in at least one call to IsAnAUGMENTINGPATH, $f$ (multiplied by some small constant) is an upper bound to the total number of computations made by LocalMM.

We state the following proposition, the proof of which appears in Section 4
Proposition 14 Let $W_{i}$ be a random variable. Let $z_{1}, z_{2}, \ldots z_{W_{i}}$ be random variables, (some possibly equal to 0 with probability 1). Assume that there exist constants $c$ and $\mu$ such that for all $1 \leq j \leq W_{i}, \operatorname{Pr}\left[z_{j} \geq \mu N\right] \leq e^{-c N}$, for all $N>0$. Then there exist constants $\mu_{i}$ and $c_{i}^{\prime}$, which depend only on $d$, such that for any $q_{i}>0$,

$$
\operatorname{Pr}\left[\sum_{j=1}^{W_{i}} z_{j} \geq \mu_{i} q_{i} \mid W_{i} \leq q_{i}\right] \leq e^{-c_{i}^{\prime} q_{i}}
$$

Using Proposition 14, we prove the following:
Proposition 15. For every $1 \leq \ell \leq 2 k-1$, there exist constants $\mu_{\ell}$ and $c_{\ell}$, which depend only on $d$ and $\epsilon$, such that for any large enough $N>0$

$$
\operatorname{Pr}\left[f_{\ell}>\mu_{\ell} N\right] \leq e^{-c_{\ell} N}
$$

Proof. The proof is by induction. For the base of the induction, we have, from Corollary 13 , that there exists an absolute constant $c_{2 k-1}$, which depends only on $d$, such that for any large enough $N>0, \operatorname{Pr}\left[X_{2 k-1}>N\right] \leq e^{-c_{2 k-1} N}$. Assume that the proposition holds for $\ell=2 k-1,2 k-3, \ldots \ell$, and we show that it holds for $\ell-2$.

Let $b_{\ell}=(\ell+2 d-2) \cdot \ell(d-1)^{\ell-1}$. From Proposition 11 we have that each call to IsAnAugmenting Path $(\ell)$ generates at most $b_{\ell}$ calls to IsPathinMIS $(\ell-2)$, and hence $b_{\ell} \cdot X_{\ell-2}$ calls to IsAnAUGMENTINGPATH $(\ell-2)$. From Corollary 13, we have that there exists an absolute constant $c$, which depends only on $d$, such that for any large enough $N>0$,

$$
\operatorname{Pr}\left[X_{\ell-2}>N\right] \leq e^{-c N}
$$

Setting $W_{\ell}=b_{\ell} f_{\ell}, f_{\ell-2}=\sum_{j=1}^{W_{i}} z_{j}, q_{i}=b_{\ell} \mu_{\ell} y_{\ell}$, and $\mu_{i}=\mu_{\ell-2} / b_{\ell} \mu_{\ell}$, and letting $c_{i}^{\prime}=c_{\ell}^{\prime} / b_{\ell} \mu_{\ell}$ in Proposition 14 implies the following:

$$
\begin{equation*}
\operatorname{Pr}\left[f_{\ell-2}>\mu_{\ell-2} y_{\ell} \mid f_{\ell} \leq \mu_{\ell} y_{\ell}\right] \leq e^{-c_{\ell}^{\prime} y_{\ell}} . \tag{1}
\end{equation*}
$$

[^4]We have

$$
\begin{align*}
\operatorname{Pr}\left[f_{\ell-2}>\mu_{\ell-2} N\right]= & \operatorname{Pr}\left[f_{\ell-2}>\mu_{\ell-2} N \mid f_{\ell} \leq \mu_{\ell} N\right] \cdot \operatorname{Pr}\left[f_{\ell} \leq \mu_{\ell} N\right] \\
& +\operatorname{Pr}\left[f_{\ell-2}>\mu_{\ell-2} N \mid f_{\ell}>\mu_{\ell} N\right] \cdot \operatorname{Pr}\left[f_{\ell}>\mu_{\ell} N\right] \\
\leq & \operatorname{Pr}\left[f_{\ell-2}>\mu_{\ell-2} N \mid f_{\ell} \leq \mu_{\ell} N\right]+\operatorname{Pr}\left[f_{\ell}>\mu_{\ell} N\right] \\
\leq & e^{-c_{\ell}^{\prime} N}+e^{-c_{\ell} N}  \tag{2}\\
= & e^{-c_{\ell-2} N},
\end{align*}
$$

where Inequality 2 stems from Inequality 1 and the induction hypothesis.
Taking a union bound over all $k$ levels immediately gives
Lemma 16. There exists a constant $c$, which depends only on $d$ and $\epsilon$, such that

$$
\operatorname{Pr}[f>c \log n] \leq 1 / n^{2}
$$

Proof (Proof of Theorem 3). Using Lemma 16, and taking a union bound over all possible queried edges gives us that with probability at least $1-1 / n$, LocalMM will require at most $O(\log n)$ queries. Therefore, for each execution of LocalMM, we require at most $O(\log n)$-independence for each conflict graph, and therefore, from Theorem 2 we require $\lceil 1 / \epsilon\rceil$ seeds of length $O\left(\log ^{3} n\right)$, which upper bounds the space required by the algorithm. The time required is upper bound by the time required to compute $r(p)$ for all the required nodes in the conflict graphs, which is $O\left(\log ^{4} n\right)$.

## 4 Combinatorial Proofs

We want to bound the total number of queries required by Algorithm LocalMM.
Let $T$ be a $d$-regular query tree. As in [2|14], we partition the interval [0,1] into $L \geq d+1$ sub-intervals: $I_{i}=$ $\left(1-\frac{i}{L+1}, 1-\frac{i-1}{L+1}\right]$, for $i=1,2, \cdots, L$ and $I_{L+1}=\left[0, \frac{1}{L+1}\right]$. We refer to interval $I_{i}$ as level $i$. A vertex $v \in T$ is said to be on level $i$ if $r(v) \in I_{i}$. Assume the worst case, that for the root of the tree, $v_{0}, r\left(v_{0}\right)=1$. The vertices on level 1 form a tree $T_{1}$ rooted at $v_{0}$. Denote the number of (sub)trees on level $i$ by $t_{i}$. The vertices on level 2 will form a forest of subtrees $\left\{T_{2}^{(1)}, \cdots, T_{2}^{\left(t_{2}\right)}\right\}$, where the total number of subtrees is at most the sum of the number of children of all the vertices in $T_{1}$. Similarly, the vertices on level $i>1$ form a forest of subtrees $F_{i}=\left\{T_{i}^{(1)}, \cdots T_{i}^{\left(t_{i}\right)}\right\}$. Note that all these subtrees $\left\{T_{i}^{(j)}\right\}$ are generated independently by the same stochastic process, as the ranks of all of the nodes in $T$ are i.i.d. random variables. Denote $f_{i}=\left|F_{i}\right|$, and let $Y_{i}=\sum_{j=1}^{i} f_{j}$. Note that $F_{i+1}$ can consist of at most $Y_{i}$ subtrees. We prove the following theorem.
Theorem 12, For any infinite query tree $T$ with bounded degree $d$, there exists a constant $c$, which depends only on $d$, such that for any large enough $N>0$,

$$
\operatorname{Pr}[|T| \geq N] \leq e^{-c N}
$$

We require the following Lemma from [14].
Lemma 17 ([14]). Let $L \geq d+1$ be a fixed integer and let $T$ be the $d$-regular infinite query tree. Then for any $1 \leq i \leq L$ and $1 \leq j \leq t_{i}$, there is an absolute constant $c$, which depends only on $d$, such that for all $N>0$,

$$
\operatorname{Pr}\left[\left|T_{i}^{(j)}\right| \geq N\right] \leq e^{-c N}
$$

We first prove the following proposition:
Proposition 18. For any infinite query tree $T$ with bounded degree d, there exist constants $\mu_{1}$ and $c_{1}$, which depend only on $d$, such that for any $1 \leq i \leq L-1$, and any $y_{i}>0$,

$$
\operatorname{Pr}\left[f_{i+1} \geq \mu_{1} y_{i} \mid Y_{i}=y_{i}\right] \leq e^{-c_{1} y_{i}}
$$

Proof. Fix $Y_{i}=y_{i}$. Let $\left\{z_{1}, z_{2}, \ldots z_{y_{i}}\right\}$ be integers such that $\forall 1 \leq i \leq y_{i}, z_{i} \geq 0$ and let $x_{i}=\sum_{i=1}^{y_{i}} z_{i}$. By Lemma 17, the probability that $F_{i+1}$ consists exactly of trees of size $\left(z_{1}, z_{2}, \ldots z_{y_{i}}\right)$ is at most $\prod_{i=1}^{y_{i}} e^{-c z_{i}}=e^{-c x_{i}}$. There are $\binom{x_{i}+y_{i}}{y_{i}}$ vectors that can realize $x_{i}{ }^{7}$ We want to bound $\operatorname{Pr}\left[f_{i+1}=\mu y_{i} \mid Y_{i}=y_{i}\right]$ for some large enough constant $\mu>0$. Letting $x_{i}=\mu y_{i}$, we bound it as follows:

$$
\begin{aligned}
\operatorname{Pr}\left[f_{i+1}=x_{i} \mid Y_{i}=y_{i}\right] & \leq\binom{ x_{i}+y_{i}}{y_{i}} e^{-x_{i}} \\
& \leq\left(\frac{e \cdot\left(x_{i}+y_{i}\right)}{y_{i}}\right)^{y_{i}} e^{-c x_{i}} \\
& =\left(\frac{e \cdot\left(\mu y_{i}+y_{i}\right)}{y_{i}}\right)^{y_{i}} e^{-c \mu y_{i}} \\
& =(e \cdot(1+\mu))^{y_{i}} e^{-c \mu y_{i}} \\
& =e^{y_{i}(-c \mu+\ln (1+\mu)+1)} \\
& \leq e^{-c^{\prime} \mu y_{i}},
\end{aligned}
$$

for some constant $c^{\prime}>0$. It follows that

$$
\begin{aligned}
\operatorname{Pr}\left[f_{i+1} \geq \mu y_{i} \mid Y_{i}=y_{i}\right] & \leq \sum_{k=\mu y_{i}}^{\infty} e^{-c^{\prime} k} \\
& \leq e^{-c_{1} y_{i}}
\end{aligned}
$$

for some constant $c_{1}>0$.
Proposition 18 immediately implies the following corollary.
Corollary 19. For any infinite query tree $T$ with bounded degree $d$, there exist constants $\mu$ and $c$, which depend only on $d$, such that for any $1 \leq i \leq L-1$, and any $y_{i}$,

$$
\operatorname{Pr}\left[f_{i+1} \geq \mu y_{i} \mid Y_{i} \leq y_{i}\right] \leq e^{-c y_{i}} .
$$

Corollary 19, which is about query trees, can be restated as follows: let $W_{i}=Y_{i}, q_{i}=y_{i}$ and $\sum_{i=1}^{W_{i}} z_{i}=f_{i+1}$. Furthermore, let $c_{i}^{\prime}=c_{1}$ and $\mu_{i}^{\prime}=\mu_{1}$ for all $i$. This notation yields the following proposition, which is unrelated to query trees, and which we used in Section 3

Proposition 14 Let $W_{i}$ be a random variable. Let $z_{1}, z_{2}, \ldots z_{W_{i}}$ be random variables (some possibly equal to 0 with probability 1). Assume that there exist constants $c$ and $\mu$ such that for all $1 \leq j \leq W_{i}, \operatorname{Pr}\left[z_{j} \geq \mu N\right] \leq e^{-c N}$, for all $N>0$. Then there exist constants $\mu_{i}$ and $c_{i}^{\prime}$, which depend only on $d$, such that for any $q_{i}>0$,

$$
\operatorname{Pr}\left[\sum_{j=1}^{W_{i}} z_{j} \geq \mu_{i} q_{i} \mid W_{i} \leq q_{i}\right] \leq e^{-c_{i}^{\prime} q_{i}} .
$$

We need one more proposition before we can prove Theorem 12 Notice that $f_{1}=\left|T_{1}\right|$.
Proposition 20. For any infinite query tree $T$ with bounded degree d, for any $1 \leq i \leq L$, there exist constants $\mu_{i}$ and $c_{i}$, which depend only on $d$, such that for and any $N>0$,

$$
\operatorname{Pr}\left[f_{i} \geq \mu_{i} N\right] \leq e^{-c_{i} N}
$$

[^5]The proof is similar to the proof of Proposition 15 We include it for completeness.
Proof. The proof is by induction on the levels $1 \leq i \leq L$, of $T$.
For the base of the induction, $i=1$, by Lemma 17, we have that there exist some constants $\mu_{1}$ and $c_{1}$ such that

$$
\operatorname{Pr}\left[f_{1} \geq \mu_{1} N\right] \leq e^{-c_{1} N}
$$

as $f_{1}=\left|T_{1}\right|$.
For the inductive step, we assume that the proposition holds for levels $1,2, \ldots, i-1$, and show that it holds for level $i$.

$$
\begin{align*}
\operatorname{Pr}\left[f_{i} \geq \mu_{i} N\right]= & \operatorname{Pr}\left[f_{i} \geq \mu_{i} N \mid Y_{i-1}<\mu_{i-1} N\right] \cdot \operatorname{Pr}\left[Y_{i-1}<\mu_{i-1} N\right] \\
& +\operatorname{Pr}\left[f_{i} \geq \mu_{i} N \mid Y_{i-1} \geq \mu_{i-1} N\right] \cdot \operatorname{Pr}\left[Y_{i-1} \geq \mu_{i-1}\right] \\
\leq & \operatorname{Pr}\left[f_{i} \geq \mu_{i} N \mid Y_{i-1}<\mu_{i-1} N\right]+\operatorname{Pr}\left[Y_{i-1} \geq \mu_{i-1}\right] \\
\leq & e^{-c N}+e^{-c_{i-1} N}  \tag{3}\\
\leq & e^{-c_{i} N}
\end{align*}
$$

for some constant $c_{i}$. Inequality 3 stems from Corollary 19 and the inductive hypothesis.
We are now ready to prove Theorem 12
Proof (Proof of Theorem[12). We would like to bound $\operatorname{Pr}\left[|T|=\sum_{i=1}^{L} f_{i} \geq \mu N\right]$. From Proposition 20, we have that for $1 \leq i \leq L$,

$$
\operatorname{Pr}\left[f_{i} \geq \mu_{i} N\right] \leq e^{-c_{i} N}
$$

A union bound on the $L$ levels gives the required result.

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## A Pseudocode for Algorithm GreedyMIS

```
Algorithm 2 - GreedyMIS - On-line MIS algorithm with input \(G=(V, E)\) and vertex permutation \(\pi\)
    \(I \leftarrow \emptyset \quad \triangleright I\) is a set of independent vertices.
    Let \(\pi=\left(v_{1}, v_{2}, \ldots, v_{n}\right)\).
    for \(i=1\) to \(n\) do
        if \(\forall u \in N\left(v_{i}\right), u \notin I\) then
            \(I \leftarrow I \cup\left\{v_{i}\right\}\).
        end if
    end for
    Output \(I\)
    \(\triangleright I\) is an MIS.
```


## B Pseudocode for Algorithm LocalMM

```
Algorithm 3-LocalMM - LCA for MM with input \(G=(V, E), e \in E\) and \(\epsilon>0\)
    Global \(\mathcal{S}=\emptyset \quad \triangleright \mathcal{S}\) is the set of seeds
    procedure \(\operatorname{Main}(G, e, \epsilon)\)
        if this is the first execution of Algorithm LocalMM then
            \((\mathcal{S}, k) \leftarrow \operatorname{Initialize}(G, \epsilon)\)
        end if
        Return \(\operatorname{IsInMAtching}(G, e, 2 k-1, \mathcal{S})\).
    end procedure
```

```
Algorithm 4 Auxiliary procedures
    procedure Initialize \((G, \epsilon) \quad \triangleright\) This is run only at the first execution
        \(k \leftarrow\lceil 1 / \epsilon\rceil\).
        for \(\ell=1,3, \ldots 2 k-1\) do
            Generate a seed \(s_{\ell}\) of length \(O\left(\log ^{3} n\right)\). \(\triangleright s_{\ell}\) is a seed for a random ordering \(\pi_{\ell}\) on all possible paths of length \(\ell\) in \(G\).
        end for
        \(\mathcal{S}=\bigcup_{\ell} \mathcal{S}_{\ell}\).
        Return \((\mathcal{S}, k)\).
    end procedure
    procedure IsInMatching \((G, e, \ell, \mathcal{S})\)
        if \(\ell=-1\) then \(\triangleright\) The empty matching
            Return false.
        end if
        \(b_{1}=\operatorname{ISINMATChing}(G, e, \ell-2, \mathcal{S})\).
        \(b_{2}=\) false.
        \(P=\{p \in G: e \in p \wedge|p|=\ell\}\)
        for all \(p \in P\) do
            if \(\operatorname{IsPathInMIS}(G, p, \ell, \mathcal{S})\) then
                \(b_{2}=\) true.
            end if
        end for
        Return \(b_{1} \oplus b_{2}\).
    end procedure
    procedure IsPathinMIS \((G, p, \ell, \mathcal{S})\)
        \(C \leftarrow \operatorname{RelevantPaths}(G, p, \ell, \mathcal{S}) . \quad \triangleright C\) is a subgraph of \(C_{M_{\ell-2}}(\ell)\)
        \(I \leftarrow\) Greedy MIS \(\left(C, \pi\left(C, s_{\ell}\right)\right)\)
        \(b=(v \in I)\)
        Return \(b\)
    end procedure
    procedure \(\operatorname{ISFreE}(G, v, \ell, \mathcal{S}) \quad \triangleright\) Checks that a vertex is free
        IsFreeVertex \(=\) true.
        for all \(u \in N(v)\) do \(\quad \triangleright\) All edges touching \(v\)
            if \(\operatorname{IsInMatching}(G,(u, v), \ell-2, \mathcal{S})\) then
                    IsFreeVertex \(=\) false.
            end if
        end for
        Return IsFreeVertex.
    end procedure
```

```
Algorithm 5 More auxiliary procedures
    procedure RelevantPaths \((G, p, \ell, \mathcal{S})\)
        Initialize \(C=\left(V_{C}, E_{C}\right) \leftarrow(\emptyset, \emptyset)\).
        if IsAnAugmentingPath \((G, p, \ell, \mathcal{S})\) then
            \(V_{C}=\{p\}\).
        else
            Return \(C\).
        end if
        while \(\exists p \in V_{C}:\left(p, p^{\prime}\right) \in E_{C}, r_{\ell}\left(p^{\prime}, s_{\ell}\right)<r_{\ell}\left(p, s_{\ell}\right)\) do
            if IsAnAugmenting Path \(\left(G, p^{\prime}, \ell, \mathcal{S}\right)\) then
                \(V_{C} \leftarrow p^{\prime}\)
                    for all \(p^{\prime \prime} \in N\left(p^{\prime}\right)\) do \(\quad \triangleright\) Edges between \(p^{\prime}\) and vertices in \(V_{C}\)
                    if \(p^{\prime \prime} \in V_{C}\) then
                    \(E_{C} \leftarrow\left(p^{\prime}, p^{\prime \prime}\right)\).
                        end if
                    end for
            end if
        end while
        Return \(C\).
    end procedure
    procedure IsAnAugmenting \(\operatorname{Path}(G, p, \ell, \mathcal{S})\)
        \(\triangleright\) Checks that \(p\) is an augmenting path.
        If \(\ell=1\) return TRUE.
                                    \(\triangleright\) all edges are augmenting paths of the empty matching
        Let \(p=\left(e_{1}, e_{2}, \ldots, e_{\ell}\right)\), with end vertices \(v_{1}, v_{\ell+1}\).
        IsPath \(=\) true.
        for \(i=1\) to \(\ell\) do
            if \(i(\bmod 2)=0\) then \(\quad \triangleright\) All even numbered edges should be in the matching
                    if \(\neg\) IsInMatching \(\left(G, e_{i}, \ell-2, \mathcal{S}\right)\) then
                        IsPath \(=\) false.
                    end if
            end if
            if \(i(\bmod 2)=1\) then \(\quad \triangleright\) No odd numbered edges should be in the matching
                if \(\operatorname{IsInMAtching}\left(G, e_{i}, \ell-2, \mathcal{S}\right)\) then
                        IsPath \(=\) false.
                    end if
            end if
        end for
        if \(\left(\neg \operatorname{ISFREE}\left(G, v_{1}, \ell, \mathcal{S}\right) \vee\left(\neg \operatorname{ISFREE}\left(G, v_{\ell+1}, \ell, \mathcal{S}\right)\right.\right.\) then
            IsPath \(=\) false. \(\quad \triangleright\) The vertices at the end should be free
        end if
        Return IsPath.
    end procedure
```


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    ${ }^{1}$ A perfect matching includes all the nodes of a bipartite graph.

[^1]:    ${ }^{2}$ Our model differs slightly from the model of [20] in that their model requires that the LCA always obeys the time and space bounds, and returns an error with some probability. It is easy to see that any algorithm which conforms to our model can be modified to conform to the model of [20] by forcing it to return an error if the time or space bound is violated.

[^2]:    ${ }^{3}$ This approach was first used by Hopcroft and Karp in [9]; however, they only applied it efficiently in the bipartite setting.
    ${ }^{4}$ Notice that the nodes of the conflict graph represent paths in $G$. Although it should be clear from the context, in order to minimize confusion, we refer to a vertex in $G$ by vertex, and to a vertex in the conflict graph by node.

[^3]:    ${ }^{5}$ Alternately, we sometimes view $r$ as a function $r:(v, s) \rightarrow[0,1]$ : Let $r^{\prime}$ be a function $r^{\prime}:(v, s) \rightarrow\left[c n^{4}\right]$, and let $f:\left[c n^{4}\right] \rightarrow$ $[0,1]$ be a function that maps each $x \in\left[c n^{4}\right]-\{1\}$ uniformly at random to the interval $\left((x-1) / c n^{4}, x / c n^{4}\right]$, and $f$ maps 1 uniformly at random to the interval $\left[0,1 / c n^{4}\right]$. Then set $r(v, s)=f\left(r^{\prime}(v, s)\right)$.

[^4]:    ${ }^{6}$ For all even $\ell$, let $f_{\ell}=0$.

[^5]:    ${ }^{7}$ This can be thought of as $y_{i}$ separators of $x_{i}$ elements.

